

Server-side Verification of Client Behavior in Cryptographic Protocols

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Abstract

Numerous exploits of client-server protocols and applications involve modifying clients to behave in ways that untampered clients would not, such as crafting malicious packets. In this paper, we demonstrate practical verification of a cryptographic protocol client’s messaging behavior as being consistent with the client program it is believed to be running. Moreover, we accomplish this without modifying the client in any way, and without knowing all of the client-side inputs driving its behavior. Our toolchain for verifying a client’s messages explores multiple candidate execution paths in the client concurrently, an innovation that we show is both specifically useful for cryptographic protocol clients and more generally useful for client applications of other types, as well. In addition, our toolchain includes a novel approach to symbolically executing the client software in multiple passes that defers expensive functions until their inputs can be inferred and concretized. We demonstrate client verification on OpenSSL to show that, e.g., Heartbleed exploits can be detected without Heartbleed-specific filtering and within seconds of the first malicious packet, and that verification of legitimate clients can keep pace with, e.g., Gmail workloads.

1 INTRODUCTION

Tampering with clients in client-server protocols or applications is an ingredient in numerous abuses. These abuses can involve exploits on the server directly, or manipulation of application state for which the client is authoritative. Examples of the former include at least ten vulnerabilities in the last two years for OpenSSL alone¹, including the high-profile Heartbleed [10] vulnerability, which enabled a tampered SSL client to extract contents of server memory. Examples of the latter are “invalid command” game cheats that permit the client greater powers or resources in the game [28].

The ideal defense would be to implement formally verified servers that incorporate all necessary input validation and application-specific checking. However, in practice, current production servers have codebases too

large to retrofit into a formally verified model. Moreover, even simple input validation is difficult to get perfect, despite extensive review, as exhibited by widely deployed security software such as implementations of Transport Layer Security (TLS). In 2014, critical vulnerabilities were discovered in all five major implementations of TLS [18], many of which turned out (after the fact) to be failures of input validation.

Since it is generally impossible to anticipate all such abuses, in this paper we explore a holistic approach to validating client behavior as consistent with a sanctioned client program. In this approach, a *verifier* monitors each client message as it is delivered to the server, to determine whether the sequence of messages received from the client so far is consistent with the program the client is believed to be running and the messages that the server has sent to the client (Fig. 1). Performing this verification is challenging primarily because inputs or nondeterministic events at the client may be unknown to the verifier, and thus, the verifier must solve for whether there exist inputs that could have driven the client software to send the messages it did. Furthermore, some of those inputs may be protected by cryptographic guarantees (private keys in asymmetric cryptography), and maliciously crafted fields may themselves be hidden by encryption, as is the case with Heartbleed.

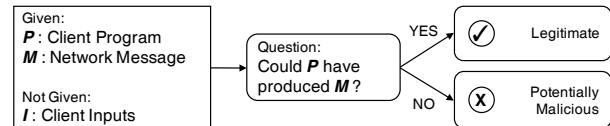


Figure 1: Abstracted client verification problem.

Our central contribution is to show that legitimate cryptographic client behavior can in fact be verified, not against a simplified protocol model but against the client program. Intuitively, we limit an attacker to only behaviors that could be effected by a legitimate client. We believe this advance to be important: in showing that messages from a client can be quickly verified as legitimate or potentially malicious, we narrow the time between zero-day exploit and detection/countermeasure to minutes or even seconds. This is significant, since in the case of Heartbleed, for example, the bug was introduced in March 2012 and disclosed in April 2014, a window of vulnerability of over two years. In fact, with latency-

¹CVE-2014-0160, CVE-2014-3512, CVE-2014-3567, CVE-2014-3513, CVE-2015-0205, CVE-2015-1787, CVE-2015-0293, CVE-2015-0292, CVE-2015-0208, CVE-2015-0291.

tolerant applications (e.g., SMTP), our defense could be installed inline as a *prevention* system rather than simply a detection system. Moreover, our technique accomplishes this verification with no vulnerability-specific configuration and, indeed, would discover client exploit attempts even prior to a vulnerability’s disclosure. It also requires no changes to client or server software.²

Following several other works in verification of client messages when some client-side values are unknown (see Sec. 2 for a discussion of prior research), the basic strategy we take is to use symbolic execution [4] to trace the client execution based on the messages received so far from the client (and the messages the server has sent to it). When the verifier, in tracing client execution, operates on a value that it does not know (a “symbolic” value), it considers all possibilities for that value (e.g., branching in both directions if a branch statement involves a symbolic variable) and records constraints on those symbolic values implied by the execution path taken. Upon an execution path reaching a message send point in the client software, the verifier reconciles the accumulated constraints on that execution path with the next message received from the client. If the path does not contradict the message, then the message is confirmed as being consistent with some valid client execution.

Our technical innovations within this direction of research are twofold.

1. Prior research on this form of client verification has primarily focused on carefully prioritizing candidate paths through the client in the hopes of finding one quickly to validate the message trace observed so far. This prioritization can itself be somewhat expensive (e.g., involving edit-distance computations on execution paths) and prone to error, in which case the verifier’s search costs grow dramatically (e.g., [7]). Here we instead use parallelism to explore candidate paths concurrently, in lieu of sophisticated path prediction or to overcome the impact of poor predictions. In Sec. 4, we describe the architecture of this parallel client-verification tool and demonstrate its ability to overcome limitations in path prediction accuracy on two online games studied in previous research. We also highlight one aspect of cryptographic protocols for which efficiently validating client-side behavior depends on being able to explore multiple execution fragments in parallel, namely execution fragments reflecting plaintexts of different sizes, when the true plaintext size is hidden by message padding (as in SSH and draft TLS 1.3). In this case, predicting the plaintext length is not possible from the ciphertext

length, by design, and so exploring different candidate lengths in parallel yields substantial savings.

2. When verifying the behavior of a client in a cryptographic protocol such as Transport Layer Security (TLS), the search for a client execution path to explain the next client message can be stymied by paths that contain cryptographic functions for which some inputs are unknown (i.e., symbolic). The symbolic execution of, e.g., the AES block cipher with an unknown message or a modular exponentiation with an unknown exponent is simply too costly. Every message-dependent branch in the AES code or modular exponentiation routine would need to be explored, and the resulting formula would be unmanageably complex. In Sec. 5 we therefore describe a multi-pass algorithm for exploring such paths, whereby user-specified “prohibitive” functions are bypassed temporarily until their inputs can be deduced through reconciliation with the client message; only then is the function explored (concretely). In cases where those inputs can never be inferred—as would be the case for an ephemeral Diffie-Hellman key, for example—the system outputs the assumption required for the verification of the client message to be correct, which can be discharged from a small whitelist of assumptions. Aside from these assumptions, our verification is sound.

We stress that our verification technique, while not completely turnkey, does not require detailed knowledge of the protocol or application being verified. For example, the specification of prohibitive functions and a matching whitelist of permissible assumptions is straightforward in our examples: the forbidden functions are simply the AES block cipher, hash functions, and elliptic curve group operations; and the whitelist includes the assumptions that a group element can be produced by raising the group generator to some power and that there exists an input that would induce the hash function to produce a given value (which are both reasonable). Aside from specifying the forbidden functions and the whitelist, the other (optional) step is to “stub out” calls to software that are irrelevant to the analysis (e.g., `printf`). For each of the contributions above, we detail the effort required by the user to utilize them.

We document the efficacy of our client verification technique by showing that it can keep pace with client messages in an interactive Gmail session running over TLS 1.2 connections. Our verification for each client-to-server TLS record takes an average of 126ms on a 3.2GHz processor. Taking into account the bursts of network activity in Gmail traffic, and given that a message cannot begin verification until all previous messages are verified, the verification of any client-to-server mes-

²In the case of validating OpenSSL client behavior, we do use a common diagnostic feature on servers: logging session keys to enable analysis of network captures.

sage completes a maximum of 14s after the time the packet was delivered over the network. We also show that our technique keeps pace with TLS connections that use message padding, a draft TLS 1.3 [9] feature that introduces costs that our parallel approach is able to overcome.

2 RELATED WORK

The most closely related work is due to Bethea et al. [1] and Cochran and Reiter [7]. These works develop algorithms to verify the behavior of (non-cryptographic) client applications in client-server settings, as we do here. Bethea et al. adopted a wholly offline strategy, owing to the expense of their techniques. Cochran and Reiter improved the method by which a verifier searches for a path through the client program that is consistent with the messages seen by the verifier so far. By leveraging a training phase and using observed messages to provide hints as to the client program paths that likely produced those messages, their technique achieved improved verification latencies but still fell far short of being able to keep pace with, e.g., highly interactive games. Their approach would not work for cryptographic protocols such as those we consider here, since without substantial protocol-specific tuning, the cryptographic protections would obscure information in messages on which their technique depends for generating these hints. In some sense, our work in Sec. 5 can be viewed as providing a method for iteratively stripping away these obfuscating effects of cryptographic fields, with the bare minimum of manual configuration or protocol-specific knowledge, and our work in Sec. 4 then dampens the impact of generated hints that are ultimately inaccurate.

Several other works have sought to verify the behavior of clients in client-server protocols. Most permit false rejections or acceptances since they verify client behavior against an abstract (and so imprecise) model of the client program (e.g., [12, 14]), versus an actual client program as we do here. Others seek exact results as we do, but accomplish this by modifying the client to send all inputs it processes to the verifier, allowing the verifier to simply replay the client on those inputs [26]. In our work, we verify completely unchanged clients and introduce no additional messaging overhead. Proxies for inferring web-form parameter constraints when a web form is served to a client, to detect parameter-tampering attacks when the form values are returned [24], also provide exact detection. However, this work addresses only stateless clients and does so without attention to cryptographically protected traffic. Our work permits stateful clients and specifically innovates to overcome challenges associated with cryptographic protocols.

Also related to our goals are numerous works focused on verifying the correctness of computations outsourced

to an untrusted cloud. Recent works in this area, surveyed by Walfish and Blumberg [27], employ advances in probabilistically checkable proofs (e.g., [15]) and/or interactive proofs (e.g., [13]) to permit a verifier to confirm (probabilistically) that an untrusted, remote party performed the outsourced computation correctly, at a cost to the verifier that is smaller than it performing the outsourced computation itself. Since we approach the problem from the opposite viewpoint of a well-resourced verifier (e.g., running with the server in a large cloud that the server owner trusts), our techniques do not offer this last property. However, ours requires no changes to the party being verified (in our case, the client), whereas these other works increase the computational cost for the party being verified (in their case, the cloud) by orders of magnitude (e.g., see [27, Fig. 5]). Another area of focus in this domain has been reducing the privacy ramifications of the additional information sent to the verifier to enable verification (e.g., [21]). Since our technique does not require changes to the messaging behavior of the application at all, our technique does not suffer from such drawbacks.

More distantly related to our work is recent progress on reducing the security of reference implementations of cryptographic protocols to underlying cryptographic assumptions (e.g., miTLS, a reference implementation of TLS in F# [2]). Whereas such works prove specified properties of client and server implementations, our work instead seeks to prove a property of the messages sent in an interaction, the property being that these messages are consistent with a specified client implementation. As such, our techniques show nothing about the intrinsic security of the client (or server) implementation itself; nevertheless, they are helpful in detecting a broad range of common exploit types in client-server protocols, as we show here. Our techniques also have the feature of being immediately deployable to existing production protocol implementations.

3 BACKGROUND AND GOALS

A client-server protocol generates messages msg_0, msg_1, \dots , some from the client and some sent by the server. Our goal is to construct a *verifier* to validate the client behavior as represented in the message sequence; the server is trusted. We assume that the client is single-threaded and that the message order reflects the order in which the client sent or received those messages, though neither of these assumptions is fundamental. Our technique is not dependent on a particular location for the verifier, though for the purposes of this paper, we assume it is near the server, acting as a passive network tap.³

³The verifier can optimistically assume that the order in which it observes the messages is the order in which the client sent or received them, and this assumption will often suffice to validate a legitimate

Borrowing terminology from prior work [7], the task of the verifier is to determine whether there exists an *execution prefix* of the client that is *consistent* with the messages msg_0, msg_1, \dots . Specifically, an execution prefix Π is a sequence of client instructions that begins at the client entry point and follows valid branching behavior in the client program. We define Π_n to be *consistent* with $msg_0, msg_1, \dots, msg_n$, if the network SEND and RECV instructions⁴ in Π_n number $n + 1$ and these network instructions match $msg_0, msg_1, \dots, msg_n$ by direction—i.e., if msg_i is a client-to-server message (respectively, server-to-client message), then the i -th network I/O instruction is a SEND (respectively, RECV)—and if the branches taken in Π_n were possible under the assumption that $msg_0, msg_1, \dots, msg_n$ were the messages sent and received. Consistency of Π_n with $msg_0, msg_1, \dots, msg_n$ requires that the conjunction of all symbolic postconditions at SEND instructions along Π_n be satisfiable, once concretized using contents of messages $msg_0, msg_1, \dots, msg_n$ sent and received on that path.

The verifier attempts to validate the sequence msg_0, msg_1, \dots incrementally, i.e., by verifying the sequence $msg_0, msg_1, \dots, msg_n$ starting from an execution prefix Π_{n-1} found to be consistent with $msg_0, msg_1, \dots, msg_{n-1}$, and appending to it an *execution fragment* that yields an execution prefix Π_n consistent with $msg_0, msg_1, \dots, msg_n$. Specifically, an *execution fragment* is a nonempty sequence of client instructions (i) beginning at the client entry point, a SEND, or a RECV in the client software, (ii) ending at a SEND or RECV, and (iii) having no intervening SEND or RECV instructions. If there is no execution fragment that can be appended to Π_{n-1} to produce a Π_n consistent with $msg_0, msg_1, \dots, msg_n$, then the search resumes by *backtracking* to find another execution prefix $\hat{\Pi}_{n-1}$ consistent with $msg_0, msg_1, \dots, msg_{n-1}$, from which the search resumes for an execution fragment to extend it to yield a $\hat{\Pi}_n$ consistent with $msg_0, msg_1, \dots, msg_n$. Only after all such attempts fail can the client behavior be declared invalid, which may take substantial time.

Determining if a program can output a given value is only semidecidable (recursively enumerable); i.e., while valid client behavior can be declared as such in finite time, invalid behavior cannot, in general. Thus, an “invalid” declaration usually comes by timeout on the verification process.⁵ However, our primary concern in this paper is verifying the behavior of *valid* clients quickly.

client even if not strictly true, particularly when the client-server protocol operates in each direction independently (as in TLS). In other cases, the verifier could in theory explore other orders when verification with the observed order fails.

⁴We abbreviate calls to POSIX `send()` and `recv()` system calls (or their functional equivalents) with the labels SEND and RECV.

⁵Nevertheless, our tool declares our tested exploit traces as invalid within several seconds; see Sec. 7.1.

4 PARALLEL CLIENT VERIFICATION

As discussed in the previous section, upon receipt of message msg_n , the verifier attempts to find an execution fragment with which to extend execution prefix Π_{n-1} (consistent with msg_0, \dots, msg_{n-1}) to create an execution prefix Π_n that is consistent with msg_0, \dots, msg_n . Doing so at a pace that keeps up with highly interactive applications remains a challenge (e.g., [7]). We observe, however, that multiple execution fragments can be explored concurrently. This permits multiple worker threads to symbolically execute execution fragments simultaneously, while coordinating their activities through data structures to ensure that they continue to examine new fragments in priority order. In this section we detail the design of our tool to do so.

While concurrent exploration of execution fragments can improve the performance of verification in any client-server application (as we show in App. A), there are specific needs for this capability for verifying the cryptographic protocols of primary interest in this paper. For example, symbolic execution tools such as the KLEE tool on which we build [6], while being designed to work with program variables whose *values* are unknown (symbolic), nevertheless require the *sizes* of those variables to be fixed. Upon observing a message msg_i that is encrypted, however, it may not be possible to determine the size of the plaintext if padding is added to the plaintext before encrypting (as in SSH and TLS 1.3)—and in some cases, this might be exactly the reason that padding was introduced (as in the case of TLS 1.3). The ciphertext length does, however, provide an upper bound on the plaintext length, and so verification can proceed by considering each possible plaintext length (up to the ciphertext length) individually. Doing so sequentially would result in a many-fold increase in verification cost, however. Instead, by considering many plaintext lengths in parallel, the increase in verification cost due to this ambiguity can be substantially dampened.

4.1 Algorithm overview

We first define the data structures used by the algorithm. A state σ represents a snapshot of execution in the symbolic virtual machine, including all constraints (path conditions) and memory objects, which includes the contents (symbolic or concrete) of registers, the stack and the heap. We use $\sigma.\text{cons}$ to represent the constraints accumulated during the execution to reach σ , and $\sigma.\text{next}$ to represent the next instruction to be executed from σ . The verifier produces state σ_n by symbolically executing the instruction sequence represented by execution prefix Π_n .

The algorithm builds and maintains a binary tree consisting of Node objects. Each node `nd` has a field `nd.path` to record a path of instructions in the client; a field `nd.state` that holds a symbolic state; children fields

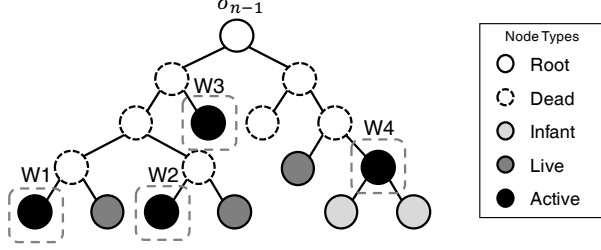


Figure 2: Example node tree.

$nd.child_0$ and $nd.child_1$ that point to children nodes, and a field $nd.saved$ that will be described in Sec. 5. The tree of nodes is rooted with a node nd holding the state $nd.state = \sigma_{n-1}$ and $nd.path = \Pi_{n-1}$. The two children of a node nd in the tree extend $nd.path$ through the next symbolic branch (i.e., branch instruction with a symbolic condition). One child node holds a state with a constraint that maintains that the branch condition implies false, and the other child node’s state holds a constraint that indicates that the branch condition is true. The algorithm succeeds by finding a fragment with which to extend Π_{n-1} to yield Π_n if, upon extending a path, it encounters a network I/O instruction that yields a state with constraints that do not contradict msg_n being the network I/O instruction’s message.

The driving goal of our algorithm is to enable concurrent exploration of multiple states in the node tree. To this end, our parallel verification algorithm uses multiple threads; it uses a single thread to manage the node tree and several worker threads, each assigned to a single node in the node tree at a time. Fig. 2 shows an example assignment of four workers to multiple nodes in a node tree. In our design and experiments, the number of worker threads $NumWorkers$ is a fixed parameter provided to the verifier. Because the verification task is largely CPU-bound, in our experience it is not beneficial to use more worker threads than the number of logical CPU cores, and in some cases, fewer worker threads than cores are necessary.

4.2 Detailed algorithm description

We will express our algorithm using standard multi-threading primitives. The keyword **spawn** indicates the creation of a child thread that will execute a named procedure until completion, at which point the child thread will terminate. The keyword **sync** denotes that the parent procedure will not proceed to the next statement until all spawned child threads have finished execution. In addition, statements involving accesses to shared data structures will be performed atomically (i.e., in a critical section), though for readability we do not include explicit designation of the critical section boundaries in our pseudocode. Also for readability, our pseudocode does not

represent backtracking.

```

100 procedure ParallelVerify( $\Pi_{n-1}, \sigma_{n-1}, msg_n$ )
101   Root  $\leftarrow$  makeNode( $\Pi_{n-1}, \sigma_{n-1}, \mathbf{true}$ )
102    $Q_R \leftarrow$  makeNodeQueue()  $\triangleright$  Nodes ready to execute
103    $Q_A \leftarrow$  makeNodeQueue()  $\triangleright$  Nodes added by workers
104   Done  $\leftarrow$  false  $\triangleright$  Instructs all threads to halt
105   Rslt  $\leftarrow$  makeNode( $\perp, \perp, \perp$ )  $\triangleright$  Filled in on success
106   spawn NodeScheduler(Root,  $Q_R, Q_A, \text{Done}$ )
107   for 1 to NumWorkers do
108     spawn VfyMsg( $msg_n, \text{Root}, Q_R, Q_A, \text{Done}, \text{Rslt}$ )
109   sync  $\triangleright$  Wait for all child threads
110   return Rslt  $\triangleright$  Return result

```

Figure 3: Main procedure for parallel client verification.

The algorithm for verifying a client-to-server message using thread-level parallelism is shown in Fig. 3. This algorithm, denoted **ParallelVerify**, takes as input the execution prefix Π_{n-1} consistent with msg_0, \dots, msg_{n-1} ; the symbolic state σ_{n-1} resulting from execution of Π_{n-1} from the client entry point on message trace msg_0, \dots, msg_{n-1} ; and the next message msg_n . Its output is $Rslt$, which holds the prefix Π_n and corresponding state σ_n in $Rslt.path$ and $Rslt.state$, respectively, if a prefix consistent with msg_0, \dots, msg_n is found. If the procedure returns with $Rslt.path = Rslt.state = \perp$, then this indicates that there is no execution prefix that can extend Π_{n-1} to make Π_n that is consistent with msg_0, \dots, msg_n . This will induce backtracking to search for another Π_{n-1} that is consistent with msg_0, \dots, msg_{n-1} , which the verifier will then try to extend to find a $\hat{\Pi}_n$ consistent with msg_0, \dots, msg_n .

The algorithm operates in a parent thread that spawns $NumWorkers + 1$ child threads; this includes one thread to manage scheduling of nodes for execution via the procedure **NodeScheduler** (not shown) and $NumWorkers$ worker threads to explore candidate execution fragments via the procedure **VfyMsg** (Fig. 4).

NodeScheduler manages the selection of node states to execute next and maintains the flow of nodes between worker threads. It receives as input two queues of nodes, a “ready” queue Q_R and an “added” queue Q_A . These queues are shared between the worker threads and the **NodeScheduler** thread. Worker threads pull nodes from Q_R and push new nodes onto Q_A . As there is only one scheduler thread and one or more worker threads producing and consuming nodes from the queues Q_R and Q_A , Q_R is a single-producer-multi-consumer priority queue and Q_A is a multi-producer-single-consumer queue. The goal of **NodeScheduler** is to keep Q_A empty and Q_R full. Nodes are in one of four possible states, either actively being explored inside **VfyMsg**, stored in Q_R , stored in Q_A , or being prioritized by **NodeScheduler**. A node at the front of Q_R is the highest priority node not currently being explored. The nodes in Q_A are child nodes that have been created by **VfyMsg** threads that

need to be prioritized by NodeScheduler and inserted into Q_R . NodeScheduler continues executing until the boolean Done is set to true by some VfyMsg thread.

Shown in Fig. 4, the procedure VfyMsg does the main work of client verification: stepping execution forward in the state σ of each node. In this figure, lines shaded gray will be explained in Sec. 5 and can be ignored for now (i.e., read Fig. 4 as if these lines simply do not exist). Like NodeScheduler, the procedure VfyMsg runs inside of a while loop until the value of Done is no longer equal to **false** (201). Recall that the parent procedure ParallelVerify spawns multiple instances of VfyMsg. Whenever there is a node on the queue Q_R , the condition on line 202 will be true and the procedure calls dequeue atomically. Note that even if $|Q_R| = 1$, multiple instances of VfyMsg may call dequeue in 203, but only one will return a node; the rest will retrieve undefined (\perp) from dequeue.

If nd is not undefined (204), the algorithm proceeds to execute the state nd.state and extend the associated path nd.path up to either the next network instruction (SEND or RECV) or the next symbolic branch (a branch instruction that is conditioned on a symbolic variable). The first case, stepping execution on a non-network / non-symbolic-branch instruction $\sigma.\text{nxt}$ (here denoted isNormal($\sigma.\text{nxt}$)), executes in a while loop on lines 206–208. The current instruction is appended to the path and the procedure execStep is called, which symbolically executes the next instruction in state σ . These lines are where the majority of the computation work is done by the verifier. The ability to concurrently step execution on multiple states is where the largest performance benefits of parallelization are achieved. Note that calls to execStep may invoke branch instructions, but these are non-symbolic branches.

In the second case, if the next instruction is SEND or RECV and if the constraints $\sigma.\text{cons}$ accumulated so far with the symbolic state σ do not contradict the possibility that the network I/O message $\sigma.\text{nxt.msg}$ in the next instruction $\sigma.\text{nxt}$ is msg_n (i.e., $(\sigma.\text{cons} \wedge \sigma.\text{nxt.msg} = \text{msg}_n) \not\Rightarrow \text{false}$, line 210), then the algorithm has successfully reached an execution prefix Π_n consistent with $\text{msg}_0, \dots, \text{msg}_n$. The algorithm sets the termination value (Done = **true**) and sets the return values of the parent function on lines 212–213: Rslt.path is set to the newly found execution prefix Π_n and Rslt.state is set to the state that results from executing it, conditioned on the last message being msg_n (denoted $[\text{execStep}(\sigma) \mid \sigma.\text{nxt.msg} \mapsto \text{msg}_n]$). All other threads of execution now exit because Done = **true** and the parent procedure ParallelVerify will return Rslt.

In the final case, (isSymbolicBranch($\sigma.\text{nxt}$)), the algorithm is at a symbolic branch. Thus, the branch condition contains symbolic variables and cannot be evalu-

```

200 procedure VfyMsg( $\text{msg}_n$ , Root,  $Q_R$ ,  $Q_A$ , Done, Rslt)
201   while  $\neg \text{Done}$  do
202     if  $|Q_R| > 0$  then
203        $\text{nd} \leftarrow \text{dequeue}(Q_R)$ 
204       if  $\text{nd} \neq \perp$  then
205          $\pi \leftarrow \text{nd.path} ; \sigma \leftarrow \text{nd.state}$ 
206         while isNormal( $\sigma.\text{nxt}$ ) do
207            $\pi \leftarrow \pi \parallel \langle \sigma.\text{nxt} \rangle$ 
208            $\sigma \leftarrow \text{execStep}(\sigma)$ 
209         if isNetInstr( $\sigma.\text{nxt}$ ) then
210           if  $(\sigma.\text{cons} \wedge \sigma.\text{nxt.msg} = \text{msg}_n) \not\Rightarrow \text{false}$  then
211             if  $(\sigma.\text{cons} \wedge \sigma.\text{nxt.msg} = \text{msg}_n) \equiv \text{nd.saved}$  then
212               Rslt.path  $\leftarrow \pi \parallel \langle \sigma.\text{nxt} \rangle$ 
213               Rslt.state  $\leftarrow$ 
214                  $[\text{execStep}(\sigma) \mid \sigma.\text{nxt.msg} \mapsto \text{msg}_n]$ 
215               Done  $\leftarrow$  true  $\triangleright$  Success!
216             else
217                $\text{nd} \leftarrow \text{clone}(\text{Root})$ 
218                $\text{nd.saved} \leftarrow \sigma.\text{cons} \wedge \sigma.\text{nxt.msg} = \text{msg}_n$ 
219               enqueue( $Q_A$ , nd)
220             else if isProhibitive( $\sigma.\text{nxt}$ ) then
221                $\text{nd.path} \leftarrow \pi \parallel \langle \sigma.\text{nxt} \rangle$ 
222                $\text{nd.state} \leftarrow \text{execStepProhibitive}(\sigma, \text{nd.saved})$ 
223               enqueue( $Q_A$ , nd)
224             else if isSymbolicBranch( $\sigma.\text{nxt}$ ) then
225                $\pi \leftarrow \pi \parallel \langle \sigma.\text{nxt} \rangle$ 
226                $\sigma' \leftarrow \text{clone}(\sigma)$ 
227                $\sigma' \leftarrow [\text{execStep}(\sigma') \mid \sigma'.\text{nxt.cond} \mapsto \text{false}]$ 
228               if  $\sigma'.\text{cons} \not\Rightarrow \text{false}$  then
229                  $\text{nd.child}_0 \leftarrow \text{makeNode}(\pi, \sigma', \text{nd.saved})$ 
230                 enqueue( $Q_A$ , nd.child0)
231                $\sigma \leftarrow [\text{execStep}(\sigma) \mid \sigma.\text{nxt.cond} \mapsto \text{true}]$ 
232               if  $\sigma.\text{cons} \not\Rightarrow \text{false}$  then
233                  $\text{nd.child}_1 \leftarrow \text{makeNode}(\pi, \sigma, \text{nd.saved})$ 
                enqueue( $Q_A$ , nd.child1)

```

Figure 4: VfyMsg procedure. Shaded lines will be explained in Sec. 5.

ated as true or false in isolation. Using symbolic execution, the algorithm evaluates both the true branch and the false branch by executing $\sigma.\text{nxt}$ conditioned on the condition evaluating to **false** (denoted $[\text{execStep}(\sigma') \mid \sigma'.\text{nxt.cond} \mapsto \text{false}]$ in line 226) and conditioned on the branch condition evaluating to **true** (230). In each case, the constraints of the resulting state are checked for consistency (227, 231), for example, using an SMT solver. If either state is consistent, it is atomically placed onto Q_A (229, 233).

4.3 Algorithm summary

Let us return to Fig. 2 from earlier, which depicts a node tree rooted at σ_{n-1} during the verification of msg_n . The node colored white with a solid outline represents the *root* node with state σ_{n-1} . The nodes colored white with dashed outlines, are the *dead* nodes and represent inter-

mediate states that no longer exist. A node is dead when it does not reach a success condition or exits the main `if` block of `VfyMsg` (starting on line 204) without generating any child nodes. Nodes colored black are the *active* nodes and are currently being explored by worker threads. Nodes colored dark gray are being prioritized by `NodeScheduler` and are still *live*. If there are worker threads that are ready to process a node, the highest priority live nodes are in Q_R . Nodes colored light gray are the *infant* nodes and are in Q_A . We can see that worker `W4` recently hit a symbolic branch condition and created two infant nodes which were added to Q_A . The other workers are likely executing lines 206–208.

While we are mainly concerned with cryptographic protocols, in App. A we show that our parallel algorithm can be highly effective in improving the speed of verifying clients in other distributed applications.

5 MULTIPASS CLIENT VERIFICATION

As shown in App. A, concurrent exploration of execution fragments can be highly beneficial to the speed of validating legitimate client behavior in non-cryptographic protocols. For verifying a cryptographic client, concurrent exploration of execution fragments can be similarly beneficial, as we will show in Sec. 7.3. Nevertheless, there remain challenges to verifying cryptographic clients that no reasonable amount of parallelization can overcome, since doing so would be tantamount to breaking some of the underlying cryptographic primitives themselves. In this section, we introduce a strategy for client verification that can overcome these hurdles for practical protocols such as TLS.

The most obvious challenge is encrypted messages. To make sense of these messages, the verifier will need to be given the symmetric session key under which they are encrypted. Fortunately, existing implementations of, e.g., OpenSSL servers, enable logging session keys to support analysis of network captures, and so we rely on such facilities to provide the session key to the verifier. Given this, it is theoretically straightforward to reverse the encryption on a client-to-server message mid-session—just as the server can—but that capability does surprisingly little to itself aid the verification of the client’s behavior. Indeed, state-of-the-art servers routinely fail to detect problems with the message sequence received from a client, as demonstrated by numerous such CVEs over the past two years in all major TLS implementations [18].

We therefore continue with our strategy of incrementally building an execution prefix Π in the client software as each message is received by the verifier to validate the client’s behavior. The verifier injects the logged session key into the execution prefix at the point where the key would first be generated by the client. Still, however, the number of execution fragments that need to be

explored in cryptographic protocol client implementations is far too large to overcome by concurrent exploration alone, when other inputs to cryptographic algorithms can be symbolic. Some of these (e.g., a message plaintext, once decrypted) could be injected by the verifier like the session key is, but in our experience, configuring where to inject what values would require substantially greater client-implementation-specific knowledge and bookkeeping than injecting just the session key does. This is in part due to the many layers in which cryptographic mechanisms are applied in modern protocols; e.g., in the TLS handshake, multiple messages are hashed to form the plaintext of another message, which is subsequently encrypted and authenticated. Even worse, other values, such as a client’s ephemeral Diffie-Hellman key, will never become available to a verifier (or server) and so cannot be injected into an execution prefix Π .

These observations motivate a design whereby the verifier is permitted to skip specified functions that would simply be too expensive to execute with symbolic inputs. Specifying such *prohibitive functions* need not require substantial client-implementation-specific or even protocol-specific knowledge; in our experience with TLS, for example, it suffices to specify basic cryptographic primitives such as modular exponentiation, block ciphers, and hash functions as prohibitive. Once specified as prohibitive, the function is skipped by the verifier if any of its inputs are symbolic, producing a symbolic result instead. Once reconciled with the message sequence msg_0, \dots, msg_n under consideration, however, the verifier can solve for some values that it was previously forced to keep symbolic, after which it can go back and verify function computations (concretely) it had previously skipped. Once additional passes yield no new information, the verifier outputs any unverified function computations (e.g., ones based on the client’s ephemeral Diffie-Hellman key) as assumptions on which the verification rests. Only if one of these assumptions is not true will our verifier erroneously accept this message trace. As we will see, however, these remaining assumptions for a protocol like TLS are minimal.

5.1 User configuration

As mentioned previously, our algorithm requires the specification of prohibitive functions. A prohibitive function is required to have no side effects other than altering its own parameters (or parameter buffers if passed by reference) and producing a return value; given the same inputs, it must produce the same results; and it must be possible to compute the sizes of all output buffers as a function of the sizes of the input buffers. A function should be specified as prohibitive if executing it on symbolic inputs induces a large number of symbolic states, due to branching that depends on input values. For ex-

ample, a physics engine might contain signal processing functions that should be marked prohibitive.

In our case studies, the prohibitive functions are cryptographic functions such as the AES block cipher or SHA-256. We stress, however, that the user need not know how these primitives are composed into a protocol. We illustrate this in App. B, where we show the user configuration needed for verifying the OpenSSL client, including the specification of the prohibitive functions.

Specifying prohibitive functions generalizes the normal procedure used by symbolic execution to inject symbolic inputs into the program. The user normally designates “user input” functions (such as `getchar`) as symbolic, so that each one is essentially replaced with a function that always returns a symbolic, unconstrained value of the appropriate size. The random number generators, client-side inputs (i.e., `stdin`), and functions that return the current time are typically so designated. The user configuration for prohibitive functions simply extends this mechanism so that some of these functions do not always return symbolic outputs, but return concrete outputs when their inputs are fully concrete.

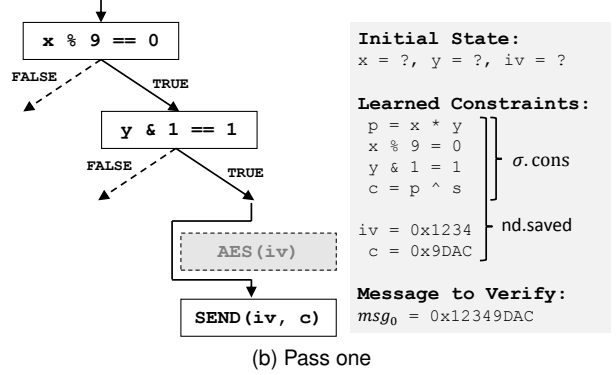
5.2 Algorithm description

The multipass verification algorithm involves changes to the `VfyMsg` procedure in Fig. 4, specifically the insertion of the shaded lines. Whenever $\sigma.\text{next}$ is a call to a prohibitive function, it is treated separately (lines 219–222), using the `execStepProhibitive` function (221). (To accomplish this, `isNormal` in line 206 now returns **false** not only for any network instruction or symbolic branch, but also for any call to a prohibitive function.) If `execStepProhibitive` receives a call $\sigma.\text{next}$ to a prohibitive function with any symbolic input buffers, it replaces the call with an operation producing fully symbolic output buffers of the appropriate size. However, if the constraints saved in `nd.saved` allow the concrete input buffer values to be inferred, then `execStepProhibitive` instead performs the call $\sigma.\text{next}$ on the now-concrete input buffers.

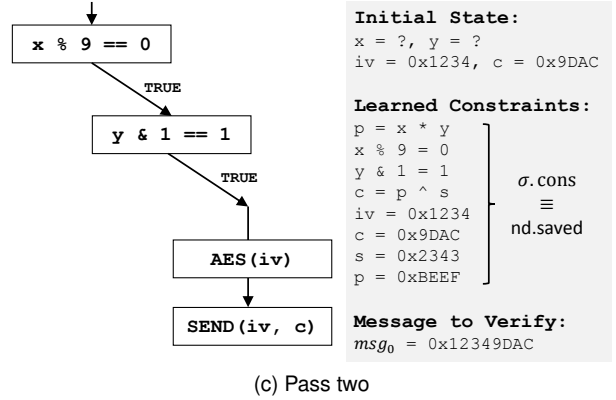
Prior to the execution path reaching a network instruction, when a call $\sigma.\text{next}$ to a prohibitive function is encountered, `nd.saved` is simply **true** as initialized (see the third argument to `makeNode` in line 101 of Fig. 3, as well as in 228 and 232), permitting no additional inferences about the values of input buffers to $\sigma.\text{next}$. After a network instruction is reached and msg_n is reconciled with the constraints $\sigma.\text{cons}$ accumulated along the path so far (210), the path constraints $\sigma.\text{cons}$ and the new constraint $\sigma.\text{next.msg} = \text{msg}_n$ are saved in `nd.saved` (217). The execution path is then replayed from the root of the binary tree (i.e., beginning from Π_{n-1} , see 216). This process repeats until an execution occurs in which nothing new is learned (i.e., $(\sigma.\text{cons} \wedge \sigma.\text{next.msg} = \text{msg}_n) \equiv \text{nd.saved}$,

```
void Client(int x, int y, int iv) {
    int p = x*y;
    if (x % 9 == 0) {
        if (y & 1 == 1) {
            int s = AES(iv);
            int c = p ^ s;
            SEND(iv, c);
        }
    }
}
```

(a) Example client code



(b) Pass one



(c) Pass two

Figure 5: Example of multipass verification on a simple client.

in 211), at which point `VfyMsg` returns as before.

5.3 Detailed walk-through

In Sec. 5.2, we concisely summarized the multipass algorithm as defined by the shaded lines in Fig. 4. We now provide a detailed walk-through of the algorithm on a segment of C code that includes both an encryption operation and a network operation (Fig. 5a), starting from entry point and following the verifier to termination.

This client multiplies two of its inputs x and y , encrypts it using a third input iv as an initialization vector, and sends both iv and the encrypted value to the server. Our tool begins with a node initialized to the client entry point and attempts to verify (by spawning one or more threads that execute `VfyMsg`) that there ex-

ist inputs x , y , and iv that would produce the output message $msg_0 = 0x12349DAC$ that was observed over the network.

The instance of `VfyMsg` that first reaches the `SEND` has, by that time, accumulated constraints $\sigma.cons$ as specified in Fig. 5b. Note, however, that it has no constraints relating s (the output of $AES(iv)$) and iv , since AES was designated as prohibitive and skipped (since iv is symbolic). After reconciling these constraints with the message $msg_0 = 0x12349DAC$, the verifier records `nd.saved`.

The verifier then re-executes from `Root` (Fig. 5c), although since it now knows $iv = 0x1234$, this time it does not skip AES . It thus computes a concrete output $s = 0x2343$ and the constraint $0x9DAC = p \wedge 0x2343$, i.e., $p = 0xBEEF$. After this second pass, the constraints in `nd.saved` are still satisfiable (e.g., $x = 0x9$, $y = 0x1537$). However, the third pass (not shown) reveals no more information, so `VfyMsg` returns the corresponding execution prefix and state at the end of the third pass.

5.4 TLS example

We illustrate the behavior of the multipass algorithm on TLS. Fig. 6 shows an abstracted subset of a TLS client implementation of AES-GCM, running on a single block of plaintext input. For clarity, the example omits details such as the implicit nonce, the server ECDH parameters, the generation of the four symmetric keys, and subsumes the tag computation into the GHASH function. But in all features shown, this walkthrough closely exemplifies the multi-pass verification of a real-world TLS client.

In Fig. 6, the outputs observed by the verifier are the client Diffie-Hellman parameter A , the initialization vector iv , the ciphertext c , and the AES-GCM tag t . The unobserved inputs are the Diffie-Hellman private exponent a , the initialization vector iv , and the plaintext p . We do assume access to the AES symmetric key k . Since the client verification is being performed on the server end of the connection, we can use server state, including the symmetric key. The verifier decides whether the observed outputs are legal, given knowledge of the program but not its inputs.

In the first pass of symbolic execution (Fig. 6a), even with knowledge of the AES symmetric key k , all prohibitive functions (ECDH, AES, GHASH) have at least one symbolic input. So, `execStepProhibitive` skips them and produces unconstrained symbolic output for each. After the first execution pass (Fig. 6b), the verifier encounters the observed client outputs. Reconciling them with the accumulated constraints $\sigma.cons$ yields concrete values for A , t , c , and iv , but not the other variables.

The verifier then begins the second pass of symbolic execution (Fig. 6c). At this point, AES and GHASH both

have concrete inputs, and therefore can be executed concretely. Note that the concrete execution of AES yields a concrete value for s , which was not previously known. At the end of the second execution pass (Fig. 6d), the verifier implicitly uses the new knowledge of s to check that there exists a p , the unobserved plaintext value, that satisfies the constraints imposed by observed output. Further passes beyond this point yield no additional information, as no further symbolic inputs to prohibitive functions can be concretized.

Note that the value of a , the client Diffie-Hellman private exponent, is never computed. The verifier thus yields as output an additional assumption that there exists an a such that $ECDH(a)$ yields values A and k . As such, we do not detect invalid curve attacks [16], for example; we discuss practical mitigations for this in Sec. 8.2.

Perhaps remarkably, no decryption mechanism is explicitly provided to the verifier. The multipass mechanism automatically recovers the plaintext for stream ciphers and counter-mode block ciphers. For other modes such as CBC, the user could provide inverse functions via an extension described in App. D.

6 IMPLEMENTATION

We have designed and implemented a prototype of our client verification technique. Our implementation is built upon a modified version of KLEE [6] and employs optimizations used in previous work [7] as well as several new performance improvements. At a high level, KLEE serves as an interpreter of LLVM assembly instructions. When an instruction has symbolic operands, the operation is stored as a symbolic expression; otherwise the operation is interpreted concretely. Note that even interpreting instructions concretely induces a significant performance penalty compared to native execution. We reduce this cost by leveraging the information provided in the user configuration of prohibitive functions to mix in native execution of calls to prohibitive functions with concrete inputs. The only additional requirement for this optimization is to provide the verifier with a native shared object that exports implementations of the prohibitive functions.

Symbolic execution of the client requires interaction with the environment through library and system calls. Our implementation inherits and extends several mechanisms from KLEE for interacting with the environment. We use a combination of the POSIX model provided by KLEE, and an expanded POSIX environment model from another system, Cloud9 [5]. We also make two key extensions to this model to support network instructions and prohibitive functions. First, during symbolic execution of the client, each `SEND` or `RECV` is intercepted and the next message in the trace under verification for a given symbolic state is processed as described

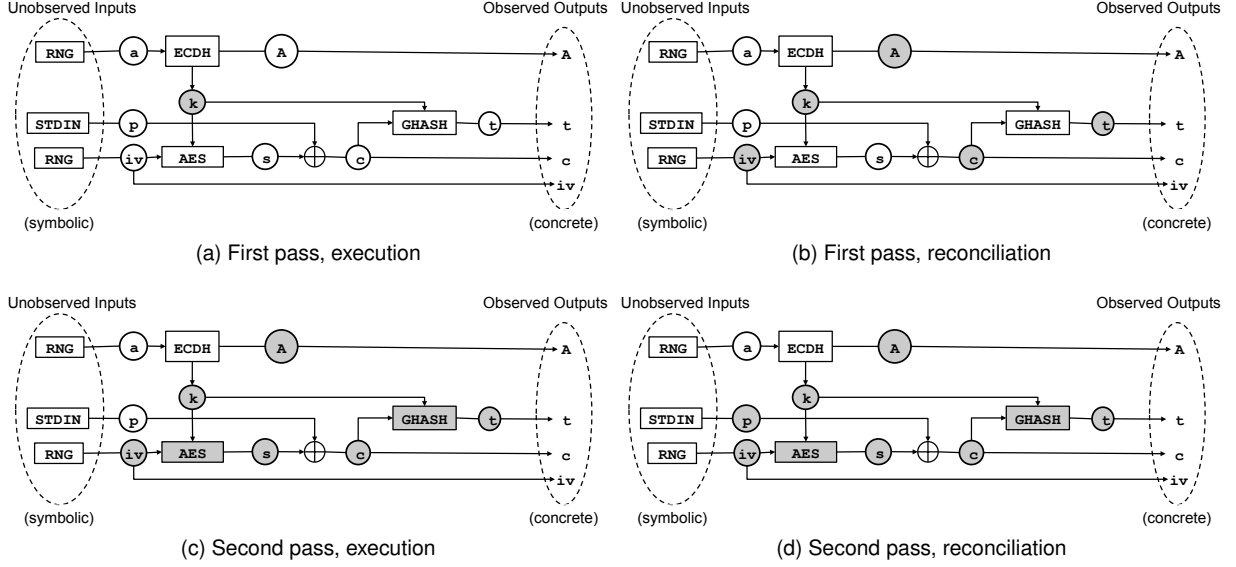


Figure 6: Multipass algorithm on a TLS client implementing an abstracted subset of AES-GCM. Rectangular blocks are prohibitive functions; circles are variables. Shaded nodes are concrete values or functions executed with concrete inputs. Unshaded nodes are symbolic values or skipped functions. In Fig. 6b and Fig. 6d, some values become concrete when $\sigma.\text{cons}$ is reconciled with $\sigma.\text{next.msg} = \text{msg}_n$ in line 210.

in Sec. 4. POSIX network calls are thus modeled to be consistent with the message trace we are verifying. Prohibitive functions are supported through an extension to the KLEE runtime model. The steps a user takes to add a prohibitive function to a given client program are straightforward. First, the user adds the function signature to a specified file in the KLEE runtime model using an API provided by the verifier. Next, the user defines the input parameters and their size, as well as the output parameters and size (as a function of the input size). The model is compiled and linked with the client program to prepare for verification. When a prohibitive function is reached during verification, the verifier uses the user-provided annotations to identify if any input memory is symbolic and, if so, the function is “skipped” and the outputs are initialized and marked as symbolic. If all inputs are concrete, then the underlying implementation is executed.

The parallel client verification algorithm (Sec. 4) requires a mechanism to concurrently execute symbolic states. We extended KLEE to support multi-threaded operation. Our changes did not significantly alter the overall architecture of KLEE, but several submodules required modification to support concurrent execution, including symbolic state “searchers”, symbolic memory management, and constraint caching and solving.

To our knowledge, no current symbolic execution engines support parallel execution in a single process via multiple threads. While other efforts have demonstrated the feasibility and performance benefits of parallelized symbolic execution engines ([5, 23, 25]), these approaches differ from ours by dividing the symbolic ex-

ecution work across multiple processes or hosts instead of across multiple threads. A multi-threaded symbolic execution engine can leverage opportunities to identify duplicate states, utilize state merging, and utilize shared constraint-solving caches. Furthermore, our architecture does not incur costs due to the latency of communication between multiple hosts; our application of symbolic execution to client verification requires high speed context switching between states. Additionally our state selection can achieve efficiencies with global knowledge of the progress of each execution path. Finally, our verifier is designed to solve SMT queries concurrently with multiple instantiations of an SMT solver, in our case STP [11].

7 EVALUATION

In this section we evaluate our implementation of the algorithms in Secs. 4–5. First, we run a single-worker verifier against two attacks on OpenSSL that represent different classes of client misbehavior. Second, we load test a single-worker verifier on a typical TLS 1.2 payload, i.e., the traffic generated by a Gmail session. Third, we increase the verification complexity to demonstrate scalability to more complex protocols with larger client state spaces, which we overcome using multiple workers. We do this by simulating verification of a TLS 1.3 draft [9] feature that permits arbitrary random padding in every packet. The OpenSSL configuration options we used are listed in App. B. All experiments were run on a system with 256GB of RAM and 3.2GHz processor cores.

Our primary measure of performance is verification *lag*. To define lag, let the verification *cost* of a message

msg_n , denoted $cost(n)$, be the wall-clock time that the verifier spends to conclude if msg_n is valid, beginning from the execution prefix Π_{n-1} produced from the successful verification of msg_0, \dots, msg_{n-1} . Since the verifier is compute-bound, $cost(n)$ is roughly the CPU time that it spends to produce Π_n from Π_{n-1} .⁶ The *completion time* for msg_n is then defined inductively as follows:

$$\begin{aligned} comp(0) &= cost(0) \\ comp(n) &= \max\{arr(n), comp(n-1)\} + cost(n) \end{aligned}$$

where $arr(n)$ is the wall-clock time when msg_n arrived at the verifier. Since the verification of msg_n cannot begin until after both (i) it is received at the verifier (at time $arr(n)$) and (ii) the previous messages msg_0, \dots, msg_{n-1} have completed verification (at time $comp(n-1)$), $comp(n)$ is calculated as the cost $cost(n)$ incurred after both (i) and (ii) are met. Finally, the *lag* of msg_n is $lag(n) = comp(n) - arr(n)$, which is our primary measure of performance.

7.1 Heartbleed and CVE-2015-0205 detection

We first evaluate our client verifier against two attacks on OpenSSL, which are meant to be illustrative of different classes of vulnerabilities that we can verify: those related to tampering with the client software to produce messages that a client could not have produced (CVE-2014-0160 Heartbleed) and a message sequence that, while correctly formatted, is impossible given a valid client state machine (CVE-2015-0205).

An OpenSSL 1.0.1f `s_server` was instantiated with standard settings, and an OpenSSL `s_client` was modified so that it would establish a TLS connection and send a single Heartbleed exploit packet. This packet had a modified length field, and when received by an OpenSSL 1.0.1f `s_server`, caused the server to disclose sensitive information from memory.

When client verification was applied to a normal client (which sends a normal Heartbeat packet), the verification lag for the (last message of the) connection was 1.7s. When it was applied to a client that sends a Heartbleed exploit, the validation process was unable to find an explanation for the contents of the packet and rejected the packet after exhausting all search paths, with a lag for the Heartbleed packet of 6.9s.

Unlike Heartbleed, CVE-2015-0205 contains only correctly formatted messages. In the certificate exchange, a good client will send a DH certificate (used to generate a pre-master secret), followed by a 0-length ClientKeyExchange message. A malicious client will send a certificate, followed by a ClientKeyExchange

message containing a DH parameter. The server will authenticate the certificate, but prefer the second message's DH parameter, allowing a malicious client to impersonate anyone whose public certificate it has obtained. We introduced this vulnerability to an OpenSSL 1.0.1d `s_server` which was instantiated with standard settings, and an OpenSSL `s_client` was modified to send a ClientKeyExchange containing a DH parameter. The server authenticated the certificate but preferred the second DH parameter.

The verification lag for the good connection was 1.3s, and the verifier rejected an attempted attack after a lag of 2.4s. This exploit illustrates the power of our technique: we not only verify whether each message is possible in isolation, but also in the context of all previous messages.

Since the tool verifies *valid* client behavior, no attack-specific configuration was required. We do not require any foreknowledge of the exploit and anticipate correct detection of other exploits requiring client tampering.

7.2 Performance evaluation: Typical TLS load

The Gmail performance test was designed to measure the lag that would result from running a single-worker verifier against typical real-world TLS traffic. The data set was a `tcpdump` capture of a three-minute Gmail session conducted in Firefox, and consisted of 21 concurrent, independent TLS sessions, totaling 3.8MB of network data. This Gmail session was performed in the context of one of the authors' email accounts and included both receiving emails and sending emails with attachments.

The verification objective of this test was to verify the TLS layer of a network connection, but not the application layer above it, such as the browser logic and Gmail web application. To simulate the client-server configuration without access to Gmail servers and private keys, we used the packet sizes and timings from the Gmail `tcpdump` to generate 21 equivalent sessions using OpenSSL `s_client` and `s_server`, such that the amount of traffic sent in each direction at any point in time matches identically with that of the original Gmail capture. Since `s_client` implements a few diagnostic features in addition to TLS (but no application layer), verifying `s_client` against these 21 sessions provides a conservative evaluation of the time required to verify the pure TLS layer.

Fig. 7 shows the bandwidth characteristics of these 21 TLS connections. As can be seen, one of the 21 TLS sessions was responsible for the vast majority of the data transferred, and almost all of the data it carried was carried from the server to the client (Fig. 7b); presumably this was a bulk-transfer connection that was involved in prefetching, attachment uploading, or other latency-insensitive tasks. The other 20 TLS sessions were utilized comparatively lightly and presumably involved in

⁶In case of backtracking—i.e., if at any point an alternate prefix $\hat{\Pi}_{n-1}$ must be produced—then the time needed to advance $\hat{\Pi}_{n-1}$ to an $\hat{\Pi}_n$ that is consistent with all of msg_0, \dots, msg_n is also accumulated into $cost(n)$.

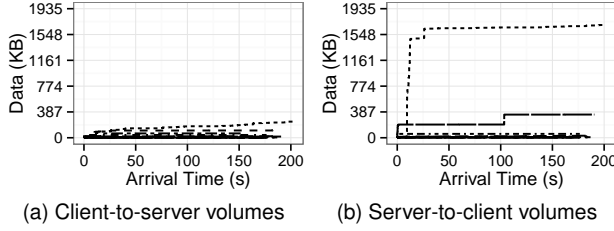


Figure 7: Cumulative data transferred for a Gmail session consisting of 21 TLS connections. Fig. 7a shows the volumes transferred in the client-to-server direction, with one line per TLS session. Fig. 7b shows the volumes transferred in the server-to-client direction.

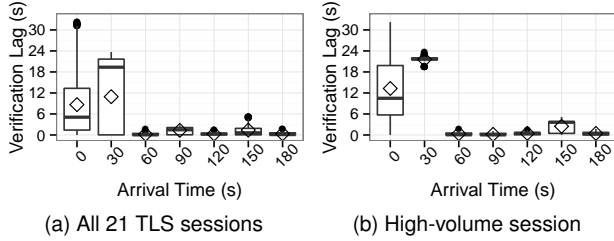


Figure 8: Verification lags in seconds for the messages in the Gmail data set. Box plot at horizontal-axis value t includes $\{lag(i) : t \leq arr(i) < t + 30s\}$. Fig. 8a shows the verification lags for all messages in all 21 TLS sessions. Fig. 8b shows the verification lags for only the messages in the high-volume session.

more latency-sensitive activities.

Fig. 8 shows the distribution of verification lag of messages, grouped by the 30-second interval in which they arrived at the verifier. Fig. 8a shows all of the messages’ verification lag, whereas Fig. 8b isolates the lag for the one high-volume flow shown in Fig. 7. It is evident from these figures that majority of the verification lag happens early on, initially up to $\sim 30s$ in the worst case. This lag coincides with an initial burst of traffic related to requests while loading the Gmail application. Another burst occurs later, roughly 160s into the trace, when an attachment was uploaded by the client. Still, even for the high-volume session (Fig. 8b), the lag is reduced to near-zero in the middle of the session and by the end of the session. The same holds true for the other 20 sessions, meaning that verification for all sessions (in parallel) can easily complete within approximately the wall-clock interval for which the sessions are active.

Fig. 9 shows verification cost as a function of mes-

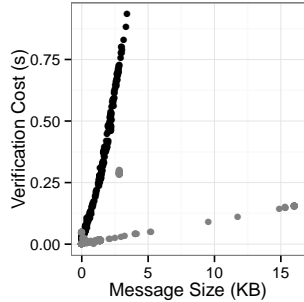


Figure 9: Message size (kilobytes) versus verification cost (seconds). Client-to-server messages (\bullet) and server-to-client handshake messages (\circ) are shown.

sage size for all 21 TLS sessions. Note that client-to-server messages tend to be costlier to verify, despite being smaller, because the verifier’s execution of the client software when processing server-to-client messages is almost entirely concrete. In contrast, the execution of the client in preparation of sending a client-to-server message tends to involve more symbolic branching. But what is most noteworthy about this plot is the linearity of the relationship between message size and verification cost, particularly for client-to-server messages. This is remarkable, given that the general problem of client verification is undecidable. In fact, it provides a simple, application-independent way to estimate the costs of verification for TLS sessions carrying payloads other than Gmail. Furthermore, assuming similar upper bounds on the sizes of messages in each direction, this predictability enables us to set a sharp timeout at which point the verifier declares the client “invalid.” Based on Fig. 9, setting the timeout at a mere 2s (verification cost) would enable the verifier to quickly detect misbehaving clients at a vanishingly small false alarm rate.

TLS-Specific Optimizations. While our goal so far in this paper has been to provide for client behavior verification with a minimum of protocol-specific tuning, a practical deployment should leverage properties of the protocol for performance. One important property of TLS (and other TCP-based protocols such as SSH) is that its client-to-server and server-to-client message streams operate independently. That is, with the exception of the initial session handshake and ending session teardown, the verifiability of client-to-server messages should be unaffected by which, if any, server-to-client messages the client has received. This gives the verifier the freedom to simply ignore server-to-client application data messages. By doing so, the verification costs for server-to-client messages, which effectively reduce to zero, do not contribute to a growing lag. The effect of this optimization on lag is shown in Fig. 10, in particular reducing the worst-case lag to around 14s. In all subsequent results, we have ignored server-to-client messages unless otherwise noted.

7.3 Stress testing: Added protocol complexity

The Gmail performance evaluation showed that verification of a typical TLS 1.2 session can be done efficiently and reliably, an advance made possible by applying a multipass methodology to cryptographic functions. In essence, once the state explosion from cryptographic functions is mitigated, the client state space becomes small enough that the verification time is primarily determined the straight-line execution speed of the KLEE symbolic interpreter. However, not all clients are guaranteed to be this simple. One good example is the draft TLS 1.3 standard [9]. In order to hide the length of the

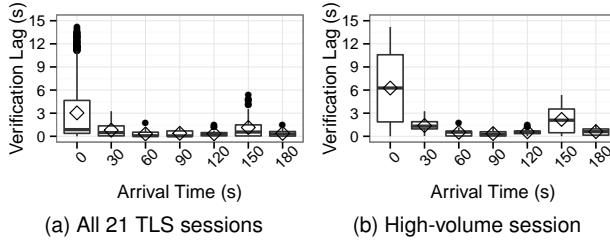


Figure 10: Verification lags in seconds for the messages in the Gmail data set, dropping server-to-client application data messages. Box plot at horizontal-axis value t includes $\{lag(i) : t \leq arr(i) < t + 30s\}$. Fig. 8a shows the verification lags for all messages in all 21 TLS sessions. Fig. 8b shows the verification lags for only the messages in the high-volume session.

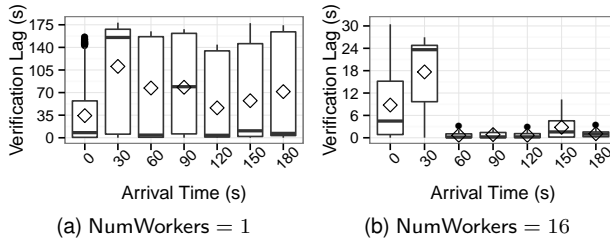


Figure 11: Verification lags in seconds for the messages in the Gmail data set when up to 128 bytes of encrypted random padding are applied, over all 21 TLS sessions. A box plot at horizontal-axis value t includes $\{lag(i) : t \leq arr(i) < t + 30s\}$. Fig. 11a shows the lag for a one-worker verifier when padding is applied. Fig. 11b shows the lag for a 16-worker verifier when padding is applied.

plaintext from an observer, implementations of TLS 1.3 are permitted (but not required) to pad an encrypted TLS record by an arbitrary size, up to maximum TLS record size. This random encrypted padding hides the size of the plaintext from any observer, whether an attacker or a verifier. In other words, given a TLS 1.3 record, the length of the input (e.g., from `stdin`) that was used to generate the TLS record could range anywhere from 0 to the record length minus header. Other less extreme examples of padding include CBC mode ciphers, and the SSH protocol, in which a small amount of padding protects the length of the password as well as channel traffic.

We thus extended our evaluation to stress test our verifier beyond typical current practice. We simulated the TLS 1.3 padding feature by modifying a TLS 1.2 client (henceforth designated as “TLS 1.2+”), so that each TLS record includes a random amount of padding up to 128 bytes⁷, applied before encryption. We then measured the performance of our verifier, ignoring server-to-client messages (except during session setup and teardown) as before.

⁷While 128 bytes of padding may seem extreme, previous work showed that an attacker could sometimes infer the website visited by encrypted HTTP connections even with substantial padding (e.g., [19]).

Figs. 11–12 show the performance of our single- and 16-worker verifiers on TLS 1.2+ with a random amount of encrypted padding. Fig. 12 shows the verification cost as a function of message size over all 21 Gmail TLS sessions and, for comparison, the cost of our baseline single-worker verifier running on a TLS 1.2 client as previously discussed in Sec. 7.2. The addition of random padding to TLS 1.2+ significantly enlarges the client state space that must be explored. When the single-worker verifier is applied to TLS 1.2+, the verification cost increases substantially compared to the TLS 1.2 baseline. The 16-worker verifier reduces the verification cost nearly back to the TLS 1.2 baseline levels. This demonstrates that the state space search is highly amenable to parallelization. Again, it is noteworthy, and perhaps surprising, that in all three cases, the verification cost is linearly related to message size.

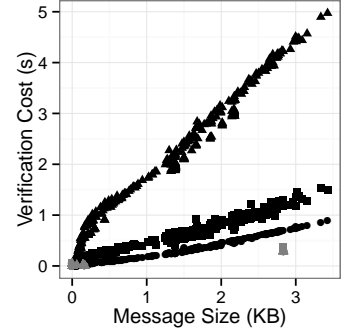


Figure 12: Message size (kilobytes) versus verification cost (seconds). Shown are the single-worker baseline verifier running on TLS 1.2 traffic (●), the single-worker verifier running on TLS 1.2+ (▲), and the 16-worker verifier running on TLS 1.2+ (■). TLS 1.2+ includes encrypted random padding. Client-to-server messages (black) and server-to-client handshake messages (gray) are also designated.

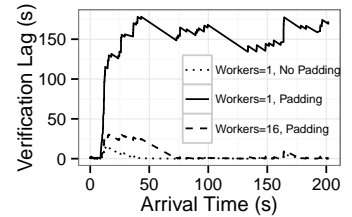


Figure 13: Verification lag in seconds for the messages in the high-volume session of the Gmail data set with ≤ 128 bytes of encrypted random padding.

Fig. 13 shows the verification lag on a TLS 1.2+ client that is sending the high-volume Gmail trace (worst case). Compared to the single-worker baseline on TLS 1.2, the single-worker verifier on TLS 1.2+ is unable to keep up with the Gmail high-volume traffic load, as its verification lag continues to increase through the end of the session at which point it is over 4 minutes behind. But with 16 workers, the initial lag (maximum of 33s lag) reduces to near-zero by the 75s mark and the verifier stays in step with the network traffic for the rest of the session. The 16-worker verifier parallelizes the search well enough to keep up with a client whose state space is two orders of magnitude larger than that of the typical TLS 1.2 client.

8 DISCUSSION

Here we discuss the applications for which our design is appropriate, and several limitations of our approach.

8.1 Applications

Suitable Application Layers. The application layer chosen for our TLS evaluation, Gmail, was a relatively strenuous one in that it exhibited both low-latency interactive behavior as well as high-volume data transfer, the latter of which was challenging as measured by verification lag. Other applications that exhibit only one behavior or the other may actually be more amenable to symbolic client verification, not only as an intrusion detection system, but as an intrusion *prevention* system.⁸ For example, the Extensible Messaging and Presence Protocol (XMPP) [22], generally used for text-based Internet messaging, is highly interactive but sends relatively small XML payloads. An average verification cost of 126ms per TLS record may be an acceptable latency for XMPP. On the other end of the spectrum are bulk data transfer applications such as the Simple Mail Transfer Protocol (SMTP). In this case, although the volume of data to be transferred can be large, the application is highly tolerant of delay. RFC 5321 recommends that the retry interval be at least 30 minutes and the give-up time be at least 4-5 days [17]. Therefore, a larger verification lag may be perfectly acceptable for TLS-protected SMTP connections.

Other Cryptographic Protocols. Perhaps due to its widespread use in various applications, TLS is one of the more complex security protocols. We believe that the client verification technique would likely generalize to other, often simpler, protocols. One example is Secure Shell (SSH), which includes both authentication and transport layer protocols [29, 30]. Indeed, SSH is likely to exhibit application behavior amenable to our verification approach. When used as a remote shell, SSH requires low latency but transfers a relatively small amount of data: key presses and terminal updates. When used for file transfer (SFTP), a large volume of data is sent, but in a mode that is relatively latency-insensitive.

8.2 Limitations

Source Code and Other TLS Implementations. Our verifier requires the client source code to generate LLVM bitcode and to designate prohibitive functions. We also

⁸Because of the need for the verifier to obtain the master secret from the server after the server receives the second client-to-server message in the TLS handshake, the verifier would need to substitute a fake master secret in place of the real one in order to verify the first two client-to-server messages, if used as an intrusion prevention system. The first *use* of the real master secret, which is not evidenced until several messages later, would then cause the verifier to backtrack, at which point the real master secret could be inserted.

require knowledge of the client configuration, such as command line parameters controlling the menu of possible cipher suites. While this work provides verification for OpenSSL-based clients, there are several other popular implementations of TLS, including Mozilla NSS, GnuTLS, Apple Secure Transport, and Microsoft Schannel. In principle, one could verify a significant portion of TLS traffic by verifying observed traffic against these five major TLS libraries, considering a client valid if its behavior is consistent with any one of them.

Client Versions. A verifier that is validating a client’s behavior against the wrong version of the client’s implementation could reject a legitimate client falsely. In App. C, we summarize our initial steps toward verifying multiple client versions far more efficiently than instantiating a separate verifier for each possible version.

Environment Modeling. While OpenSSL `s_client` has relatively few interactions with the environment, other clients may interact with the environment extensively. For example, *XPilot* interacts with the windowing system, and SSH reads `/etc/passwd`, the `.ssh/` directory, redirects standard file descriptors, etc. The KLEE [6] and Cloud9 [5] POSIX runtimes serve as good starting points, but some environment modeling is likely to be necessary for each new type of client. This one-time procedure per client is probably unavoidable.

Manual Choice of Prohibitive Functions. We currently choose the set of prohibitive functions manually. While the choice of cryptographic hash functions, public key algorithms, and symmetric cipher primitives may be relatively obvious to a security researcher, this may not be true for a typical software developer.

Prohibitive Function Assumptions. When prohibitive functions are initially skipped but eventually executed concretely, verification soundness is preserved. If a prohibitive function is never executed concretely (e.g., due to asymmetric cryptography), this introduces an assumption; e.g., in the case of ECDH, a violation of this assumption could yield an invalid curve attack [16]. In a practical deployment, the user designating a prohibitive function should also designate predicates on the function’s output (e.g., the public key is actually a group element) that are specified by the relevant NIST or IETF standards as mandatory server-side checks [20] (which would have prevented the Jager et al. attack [16]). In our tool, these predicates could be implemented via lazy constraint generation (see App. D), or as a `klee_assume` for simple predicates. Of course, care must be taken not to turn our verifier into an oracle that enables Bleichenbacher-type attacks [3].

IDS/IPS Deployment and Denial of Service. Our current tool takes as input a recorded network log. To de-

ploy it as an intrusion detection system (IDS), it would need to be integrated with a passive network tap. For latency-tolerant applications, it could even be deployed as an intrusion prevention system (IPS), where it acts as a firewall that delivers only verified packets to the server. In both modes, but especially the latter, we must mitigate a potential denial of service (DoS) attack. To do so, we leverage the strong linear relationship between verification cost and message size in two ways. (1) Impose a hard upper bound on verifier time per packet, and declare all packets that exceed the time budget invalid. Since legitimate packets finish within a few seconds, the bound can be easily set such that the false alarm rate is negligible. (2) Given a fixed CPU time budget, precisely compute the amount of traffic that can be verified. The operator can then allocate verifiers according to the threat profile, e.g., assigning verifiers to high-priority TLS sessions or ones from networks with poor reputation (e.g., [8]). The IDS/IPS will then degrade gracefully when total traffic bandwidth exceeds the verification budget.

9 CONCLUSION

We showed that it is possible to efficiently verify that the messaging behavior of an untrusted cryptographic client is consistent with legitimate client code. Our technical contributions are twofold. First, we described a parallel client verification engine that supports concurrent exploration of paths in the client software to explain a sequence of observed messages. This innovation is both generally useful for client verification and specifically useful for verifying cryptographic clients, e.g., due to ambiguities arising from message padding hidden by encryption. Second, we developed a multipass verification strategy that enables verification of clients whose code contains cryptographic functions, which typically pose major challenges to symbolic execution. We demonstrated our verifier by showing that it detects two attacks on OpenSSL that represent two classes of client misbehavior: those that produce malformed messages (e.g., Heartbleed), and those that leverage correctly formatted messages that are nevertheless impossible to observe from a valid client. In addition, we showed that our verifier can keep pace with a typical TLS load (Gmail), running over both OpenSSL TLS 1.2 and a more complex simulation of TLS 1.3. We believe our contributions to be significant in reducing the detection time of a nontrivial class of protocol exploits from years to seconds, with no prior knowledge of the exploit.

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A DEMONSTRATION OF PARALLELIZATION ON CLIENT-SERVER GAMES

Though our focus in this paper is on cryptographic protocols, the benefits of the parallel verification architecture described in Sec. 4 extend to other protocols, as well. To demonstrate this, we employ two game clients studied in previous work [7], namely *XPilot* and *TetriNet* (specifically *XPilot NG* v4.7.2 and *TetriNet* v0.11). The *XPilot* client consists of roughly 100,000 SLOC. Beyond this, the scope of symbolic execution included all needed libraries except `Xlib`, whose functions were replaced with minimal stubs, so that the game could be run without display output. Moreover, `uClibc` was used in lieu of the GNU C library. The *TetriNet* client is 5000 SLOC. As in *XPilot*, the scope of symbolic execution also included all needed libraries, though again the display output library (`ncurses`) was disabled using minimal stub functions and `uClibc` was used in place of the GNU C library. The experiments shown here were run on a system with 256GB of RAM and 3.2GHz processor cores.

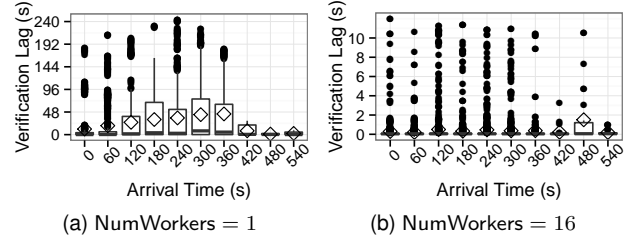


Figure 14: Verification lag of legitimate *TetriNet* traces. Box plot at horizontal-axis value t includes $\{lag(i) : t \leq arr(i) < t + 60s\}$ in 20 traces, each starting at time $t = 0$. “ \diamond ” shows the average.

Execution fragments were prioritized by *SelectNode* as described in prior work [7].

Fig. 14 shows the distribution of verification lag per message, binned into 60s bins, across 20 *TetriNet* traces. The boxplot labeled t shows the distribution of verification lags for messages that arrived between times t and $t + 60s$ in the 20 traces. In each boxplot, the “box” shows the first, second (median) and third quartiles, and its whiskers extend to cover points within ± 1.5 times the interquartile range. Additional outlier points are shown as dots. Overlaid on each boxplot is a diamond (\diamond) that shows the average of the data points. In the single worker configuration (Fig. 14a), verification lags behind message arrival times by more than 200s in the worst case. In contrast, with 16 workers (Fig. 14b), verification is able to keep pace with gameplay and never accumulates lag over the course of verification. Even if verification falls behind at some point in the game, it always catches up because of the gap between message arrival times. As such, the verifier should need only a fixed sized buffer of network messages to manage a long-running verification session.

In Fig. 15, the verification lags for *XPilot* are shown for two worker configurations across 40 message traces. Despite a mean verification cost of 75ms when using a single worker thread, the fast pace of *XPilot* makes it difficult for verification to keep pace with the game (Fig. 15a). However, by increasing the number of worker threads, we can see that in the 8-worker configuration (Fig. 15b), verification lag never significantly falls behind and could use only a fixed buffer of messages for verification in long-running sessions. In this case, there is little additional improvement gained by moving to 16 workers.

B TLS EXPERIMENTAL SETUP

In Sec. 7, we applied our client verification algorithm to OpenSSL, a widely used implementation of Transport Layer Security (TLS) with over 400,000 lines of code. We configured our OpenSSL client with one of the currently preferred cipher suites, namely `TLS_ECDHE_ECDSA_WITH_AES_128_GCM_SHA256`.

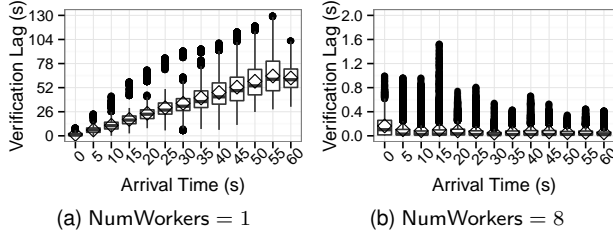


Figure 15: Verification lag of legitimate *XPilot* traces. Box plot at horizontal-axis value t includes $\{lag(i) : t \leq arr(i) < t + 5s\}$ in 40 traces, each starting at time $t = 0$. “ \diamond ” shows the average.

- Key exchange: Ephemeral Elliptic-Curve Diffie Hellman (ECDHE) signed using the Elliptic Curve Digital Signature Algorithm (ECDSA)
- Symmetric Encryption: 128-bit Advanced Encryption Standard (AES) in Galois/Counter Mode
- Pseudorandom function (PRF) built on SHA-256

Since our goal was to verify the TLS layer and not the higher-layer application, in our experiment we took advantage of the OpenSSL `s_client` test endpoint. This client establishes a fully functional TLS session, but allows arbitrary application-layer data to be sent and received via `stdin` and `stdout`, similar to the `netcat` tool. Verifying that network traffic is consistent with `s_client` is roughly equivalent to verifying the TLS layer alone.

The OpenSSL-specific user configuration for verification consisted of the following:

1. Configuring the following OpenSSL functions as prohibitive: `AES_encrypt`, `ECDH_compute_key`, `EC_POINT_point2oct`, `EC_KEY_generate_key`, `SHA1_Update`, `SHA1_Final`, `SHA256_Update`, `SHA256_Final`, `gcm_gmult_4bit`, `gcm_ghash_4bit`
2. Configuring `tls1_generate_master_secret` as the function to be replaced by server-side computation of the symmetric key.
3. (Optional) Declaring `EVP_PKEY_verify` to be a function that always returns success. Performance optimization only.

The user configuration for OpenSSL, comprising declarations of prohibitive functions and their respective input/output annotations, consisted of 138 lines of C code using our API. Fig. 16 shows an example prohibitive function declaration for the AES block cipher. In this macro, we declare the function signature, which comprises the 128-bit input buffer `in`, the 128-bit output buffer `out`, and the symmetric key data structure, `key`, which contains the expanded round keys. Both `in` and

```
DEFINE_MODEL(void, AES_encrypt,
             const unsigned char *in,
             unsigned char *out,
             const AES_KEY *key)
{
    SYMBOLIC_CHECK_AND_RETURN(
        in, 16,
        out, 16, "AESBlock");
    SYMBOLIC_CHECK_AND_RETURN(
        key, sizeof(AES_KEY),
        out, 16, "AESBlock");
    CALL_UNDERLYING(AES_encrypt,
        in, out, key);
}
```

Figure 16: Example prohibitive function declaration.

`key` are checked for symbolic data. If either buffer contains symbolic data, `out` is populated with unconstrained symbolic data, and the macro returns without executing any subsequent lines. Otherwise, the underlying (concrete) AES block cipher is called.

In a pure functional language or an ideal, strongly typed language, the prohibitive function specifications could in principle be generated automatically from the function name alone. Unfortunately, in C, the memory regions representing input and output may be accessible only through pointer dereferences and type casts. This is certainly true of OpenSSL (e.g., there is no guarantee that the `AES_KEY` struct does not contain pointers to auxiliary structs). Therefore, for each prohibitive function, the user annotation must explicitly define the data layout of the input and output.

The domain knowledge required for the first two configuration steps is minimal, namely that current TLS configurations use the above cryptographic primitives in some way, and that a symmetric key is generated in a particular function. The domain knowledge necessary for the third configuration step is that TLS typically uses public key signatures only to authenticate the server to the client, e.g., via the Web PKI. The server itself generates the signature that can be verified via PKI, and so the verifier knows that the chain of signature verifications is guaranteed to succeed. Moreover, this optimization generalizes to any protocol that uses a PKI to authenticate the server to an untrusted client.

C MULTI-VERSION VERIFIERS

When the version of the client software used by the verifier is different from the version of the client software being run by a legitimate client, it is possible for the verifier to falsely accuse the client as being invalid. This poses a challenge for verification when the client version is not immediately evident to the verifier. For example, OpenSSL does not communicate the minor version number of its client to the server, and hence our

verifier would be in the dark about this version number. The possibility for false accusations here is real: we confirmed, for example, that a verifier for OpenSSL client version 1.0.1e can fail if used to verify traffic for OpenSSL client version 1.0.1f, and vice versa. This occurs because, for example, the changes from 1.0.1e to 1.0.1f included removing MD5 from use and removing a timestamp from a client nonce, not to mention numerous platform-specific adaptations and bug fixes. In total, 1.0.1f involved changes to 102 files amounting to 1564 insertions and 997 deletions (according to `git`), implemented between Feb 11, 2013 and Jan 6, 2014.

The immediate solution to this problem is to instantiate a verifier for every possible version that a legitimate client might be using. By running these verifiers in parallel on the message trace from the client, the trace can be considered valid as long as one verifier remains accepting of the trace. Running many verifiers in parallel incurs considerable expense, however.

An alternative approach to address this issue is to create a single verifier that verifies traffic against several versions simultaneously — a *multi-version verifier* — while amortizing verification costs for their common code paths across all versions. To illustrate the potential savings, we constructed a multi-version verifier for both 1.0.1e and 1.0.1f by manually assembling a “unioned client” of these versions, say client “1.0.1ef”. In client 1.0.1ef, every difference in the code between client 1.0.1e and client 1.0.1f is preceded by a branch on version number, i.e.,

```
if (strcmp(version, "1.0.1e") == 0) {
    /* 1.0.1e code here */
} else {
    /* 1.0.1f code here */
}
```

We then provided this as the client code to the verifier, marking `version` as a symbolic variable. It is important to note that once the client messages reveal behavior that is consistent with only one of 1.0.1e and 1.0.1f, then `version` will become concrete, causing the verifier to explore only the code paths for that version; as such, the verifier still allows only “pure 1.0.1e” or “pure 1.0.1f” behavior, not a combination thereof.

As shown in Table 1, the single-worker costs (specifically, $\sum_i \text{cost}(i)$) of verifying 1.0.1e traffic with a 1.0.1ef verifier and of verifying 1.0.1f traffic with a 1.0.1ef verifier were both within 4% of the costs for verifying with a 1.0.1e and 1.0.1f verifier, respectively. (For these tests, we used the same Gmail traces used in Sec. 7.) In fact, we do not include lag graphs for the multi-version verifier here (analogous to those in Sec. 7) because they are visibly indistinguishable from

Measurement	Network Trace	
	1.0.1e	1.0.1f
Symbolic branches	32.3%	32.0%
SMT solver queries	7.8%	7.6%
Verification cost	3.0%	3.3%

Table 1: Percentage overhead incurred by verification with a “union” 1.0.1ef client instead of the matching client. Verification cost (time) is defined in Sec. 7.

those for single-version verifiers. As can be seen from Table 1, despite a 32% increase in symbolic branches that in turn drives a 7% increase in SMT solver queries, the overall cost increases very little. This implies that despite an increase in the number of code path “options” comprising the union of two versions of client code, the incorrect paths die off quickly and contribute relatively little to total verification cost.

While a demonstration of the efficiency of a multi-version verifier for only two versions of one codebase, we believe this result suggests a path forward for verifying clients of unknown versions much more efficiently than simply running a separate verifier for each possibility. We also anticipate that multi-version verifiers can be constructed automatically from commit logs to repositories, a possibility that we hope to explore in future work.

D EXTENSION: LAZY CONSTRAINT GENERATORS

There are several potentially useful extensions to our client verification algorithm that we are considering for future development. Here we highlight one, namely *lazy constraint generators* to accompany the designation of prohibitive functions. Since a function, once specified as prohibitive, will be skipped by the verifier until its inputs are inferred concretely, the verifier cannot gather constraints relating the input and output buffers of that function until the inputs can be inferred via other constraints. There are cases, however, where introducing constraints relating the input and output buffers once some subset of them are inferred concretely would be useful or, indeed, is central to eventually inferring other inputs concretely.

Perhaps the most straightforward example arises in symmetric encryption modes that require the inversion of a block cipher in order to decrypt a ciphertext (e.g., CBC mode). Upon reaching the client SEND instruction for a message, the verifier reconciles the observed client-to-server message msg_n with the constraints $\sigma.\text{cons}$ accumulated on the path to that SEND; for example, suppose this makes concrete the buffers corresponding to outputs of the encryption routine. However, because the block cipher was prohibitive and so skipped, constraints relating the input buffers to those output buffers were not recorded, and so the input buffers remain unconstrained by the (now concrete) output buffers. Moreover, a second

pass of the client execution will not add additional constraints on those input buffers, meaning they will remain unconstrained after another pass.

An extension to address this situation is to permit the user to specify a lazy constraint generator along with designating the block cipher as prohibitive. The lazy constraint generator would simply be a function from some subset of the input and output buffers for the function to constraints on other buffers. The generator is “lazy” in that it would be invoked by the verifier only after its inputs were inferred concretely by other means; once invoked, it would produce new constraints as a function of those values. In the case of the block cipher, the most natural constraint generator would be the inverse function, which takes in the key and a ciphertext and produces the corresponding plaintext to constrain the value of the input buffer.

Our OpenSSL case study in Sec. 5.4 does not require this functionality since in the encryption mode used there, the ciphertext and plaintext buffers are related by simple exclusive-or against outputs from the (still prohibitive) block cipher applied to values that can be inferred concretely from the message. So, once the inputs to the block cipher are inferred by the verifier, the block cipher outputs can be produced concretely, and the plaintext then inferred from the concrete ciphertexts by exclusive-or.